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2017

Kontinen , J , Müller , J-S , Schnoor , H & Vollmer , H 2017 , ' Modal Independence Logic ' , þÿ J o u r n a l o f L o g i c a n d C o m p u t a t i o n , v o l . 2 7 , n o . 5 , p p . 1 3 3 3 1 3 5 2 . h t

http://hdl.handle.net/10138/214807
https://doi.org/10.1093/logcom/exw019

acceptedVersion

Modal Independence Logic

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Abstract

This paper introduces modal independence logic MIL, a modal logic that can explicitly talk about independence among propositional variables. Formulas of MIL are not evaluated in worlds but in sets of worlds, so called *teams*. In this vein, MIL can be seen as a variant of Väänänen's modal dependence logic MDL. We show that MIL embeds MDL and is strictly more expressive. However, on singleton teams, MIL is shown to be not more expressive than usual modal logic, but MIL is exponentially more succinct. Making use of a new form of bisimulation, we extend these expressivity results to modal logics extended by various generalized dependence atoms. We demonstrate the expressive power of MIL by giving a specification of the anonymity requirement of the *dining cryptographers* protocol in MIL. We also study complexity issues of MIL and show that, though it is more expressive, its satisfiability and model checking problem have the same complexity as for MDL.

Keywords: dependence logic, team semantics, independence, expressivity over finite models, computational complexity.

1 Introduction

The concept of independence is ubiquitous in many scientific disciplines such as experimental physics, social choice theory, computer science, and cryptography. Dependence logic D, introduced by Jouko Väänänen in [16], is a new logical framework in which various notions of dependence and independence can be formalized and studied. Dependence logic extends first-order logic by so called dependence atoms

$$=(x_1, \ldots, x_{n-1}, x_n),$$

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expressing that the value of the variable x_n depends (only) on the values of $x_1, \ldots, x_{n-1}$, in other words, that x_n is functionally dependent of $x_1, \ldots, x_{n-1}$. Of course, such a dependency does not make sense when talking about single assignments; therefore dependence logic formulas are evaluated for so called *teams*, i. e., sets of assignments. A team can for example be a relational database table, a collection of plays of a game, or a set of agents with features. It is this *team semantics*, together with dependence atoms, that gives dependence logic its expressive power: it is known that D is as expressive as Σ^1_1, that is, the properties of finite structures that can be expressed in dependence logic are exactly the NP-properties.

Independence logic [8] is a variant of dependence logic defined in terms of independence atoms

$$(x_1, \ldots, x_\ell) \perp_{(z_1, \ldots, z_m)} (y_1, \ldots, y_n),$$

instead of dependence atoms. The intuitive meaning of the atom is that, when the values of the variables $z_1, \ldots, z_m$ are fixed, knowing the values of $x_1, \ldots, x_\ell$ does not tell us anything new about the values of $y_1, \ldots, y_n$. Independence logic has turned out to be a very interesting variant of dependence logic. In particular, independence atoms correspond to a widely studied class of database dependencies called embedded multivalued dependencies.

In a slightly later paper Väänänen [17] introduced dependence atoms into (propositional) modal logic. Here, teams are sets of worlds, and a dependence atom $=(p_1, \ldots, p_{n-1}, p_n)$ holds in a team T if there is a Boolean function that determines the value of p_n from those of $p_1, \ldots, p_{n-1}$ in all worlds in T. The so obtained modal dependence logic MDL was studied from the point of view of expressivity and complexity in [15].

In this article we introduce a modal variant of independence logic called *modal independence logic*, MIL, extending the formulas of modal logic ML by *independence* atoms

$$(p_1, \ldots, p_\ell) \perp_{(r_1, \ldots, r_m)} (q_1, \ldots, q_n),$$

the meaning of which is that the propositional sequences $\vec{p}$ and $\vec{q}$ are independent of each other for any fixed value of $\vec{r}$. In modal independence logic, dependencies between propositions can be expressed, and thus, analogously to the first-order case, MDL can be embedded as a sublogic into MIL, and it is easy to see that MIL is strictly more expressive than MDL.

The aim of this paper is to initiate a study of the expressiveness and the computational complexity of modal independence logic. For this end, we first study the computational complexity of the satisfiability and the model checking problem for MIL. We show that, though MIL is more expressive than MDL, the complexity of these decision problems stays the same, i. e., the satisfiability problem is complete for nondeterministic exponential time (NEXP-complete, [15]) and the model checking problem is NP-complete [6]. In order to settle the complexity of satisfiability for MIL, we give a translation of MIL-formulas

to existential second-order logic formulas the first-order part of which is in the Gödel–Kalmár–Schütte prefix class. Our result then follows from the classical result that the satisfiability problem for this prefix class is NEXP-complete [3]. We will also show that the same upper bound on satisfiability can be obtained for a whole range of variants of MIL via the notion of a generalized (modal) dependence atom (a notion introduced in the first-order framework in [13]).

The expressive power of MDL was first studied by Sevenster [15], where he showed that MDL is equivalent to ML on singleton teams. In this paper we prove a general result showing that MIL, and in fact any variant of it whose generalized dependence atoms are FO-definable, is bound to be equivalent to ML over singleton teams. Interestingly, it was recently shown in [5] that a so-called extended modal dependence logic EMDL is strictly more expressive than MDL even on singletons.

To demonstrate the potential applications of MIL, we consider the *dining cryptographers* protocol [4], a classic example for anonymous broadcast which is used as a benchmark protocol in model checking of security protocols [1]. We show how the anonymity requirement of the protocol can be formalized in modal independence logic, where—unlike in the usual approaches using epistemic logic—we do not need to use the Kripke model's accessibility relation to encode knowledge, but to express the "possible future" relation of branching-time models. In addition to demonstrating MIL's expressivity, we also derive a succinctness result from our modeling of the dining cryptographers: While MIL and ML are equally expressive on singletons, MIL is exponentially more succinct.

2 Modal Independence Logic

Definition 2.1 The syntax of *modal logic* ML is inductively defined by the following grammar in extended Backus–Naur form:

$$\phi ::= p \mid \overline{p} \mid \phi \wedge \phi \mid \phi \vee \phi \mid \Diamond\phi \mid \Box\phi.$$

The syntax of *modal dependence logic* MDL is defined by

$$\phi ::= p \mid \overline{p} \mid =(\vec{q}, p) \mid \phi \wedge \phi \mid \phi \vee \phi \mid \Diamond\phi \mid \Box\phi,$$

where p is a propositional variable and $\vec{q}$ a sequence of propositional variables.

The syntax of *modal independence logic* MIL is defined by

$$\phi ::= p \mid \overline{p} \mid \vec{p} \perp_{\vec{r}} \vec{q} \mid \phi \wedge \phi \mid \phi \vee \phi \mid \Diamond\phi \mid \Box\phi,$$

where p is a propositional variable and $\vec{p}, \vec{r}, \vec{q}$ are sequences of propositional variables. The sequence $\vec{r}$ may be empty.

It is worth noting that the negation of dependence logic, and that of MDL, is not the classical negation but a so-called "game theoretic" negation that still satisfies the usual De Morgan laws. As now customary with team semantics, we think of negation as a defined operation and restrict attention to formulas in

which negation only appears in front of proposition symbols (formulas $\neg{=}(\vec{q},p)$ are logically equivalent to falsum and hence can also be dispensed without loss of generality [17]). A more detailed account of the role of negation in dependence logic can be found in [12].

A Kripke structure is a tuple $\mathcal{M} = (W, R, \pi)$, where W is a non-empty set of worlds, R is a binary relation over W and $\pi\colon W \to \mathcal{P}(\mathrm{V})$ is a labeling function for a set V of propositional variables. A *team* is a (possibly empty) set $T \subseteq W$. As usual, in a Kripke structure $\mathcal{M}$ the set of all successors of $T \subseteq W$ is defined as $R(T) = \{s \in W \mid \exists s' \in T : (s', s) \in R\}$. Furthermore we define $R\langle T\rangle = \{T' \subseteq R(T) \mid \forall s \in T\ \exists s' \in T' : (s, s') \in R\}$, the set of legal successor teams.

Definition 2.2 Let $\vec{p} = (p_1, \ldots, p_n)$ be a sequence of variables and w, w' be worlds of a Kripke model $\mathcal{M} = (W, R, \pi)$. Then w and w' are equivalent under π over $\vec{p}$, denoted by $w \equiv_{\pi,\vec{p}}\, w'$, if the following holds:

$$\pi(w) \cap \{p_1, \ldots, p_n\} = \pi(w') \cap \{p_1, \ldots, p_n\}.$$

Definition 2.3 (Semantics of ML, MDL, and MIL) Let $\mathcal{M} = (W, R, \pi)$ be a Kripke structure, T be a team over $\mathcal{M}$ and ϕ be a formula. The semantic evaluation (denoted as $\mathcal{M}, T \models \phi$) is defined inductively as follows.

$$\begin{array}{lcl}
\mathcal{M}, T \models p & \Leftrightarrow & \forall w \in T\colon p \in \pi(w) \\
\mathcal{M}, T \models \overline{p} & \Leftrightarrow & \forall w \in T\colon p \notin \pi(w) \\
\mathcal{M}, T \models \phi_1 \wedge \phi_2 & \Leftrightarrow & \mathcal{M}, T \models \phi_1 \text{ and } \mathcal{M}, T \models \phi_2 \\
\mathcal{M}, T \models \phi_1 \vee \phi_2 & \Leftrightarrow & \exists T_1, T_2\colon T_1 \cup T_2 = T, \mathcal{M}, T_1 \models \phi_1 \text{ and } \mathcal{M}, T_2 \models \phi_2 \\
\mathcal{M}, T \models \Diamond\phi & \Leftrightarrow & \exists T' \in R\langle T\rangle\colon \mathcal{M}, T' \models \phi \\
\mathcal{M}, T \models \Box\phi & \Leftrightarrow & \mathcal{M}, R(T) \models \phi \\
\mathcal{M}, T \models {=}(\vec{q}, p) & \Leftrightarrow & \forall w, w' \in T\colon w \equiv_{\pi,\vec{q}}\, w' \text{ implies } w \equiv_{\pi,p} w' \\
\mathcal{M}, T \models \vec{p}_1 \perp_{\vec{q}} \vec{p}_2 & \Leftrightarrow & \forall w, w' \in T\colon w \equiv_{\pi,\vec{q}}\, w' \text{ implies } \exists w'' \in T\colon \\
 & & w'' \equiv_{\pi,\vec{p}_1} w \text{ and } w'' \equiv_{\pi,\vec{p}_2} w' \text{ and } w'' \equiv_{\pi,\vec{q}}\, w
\end{array}$$

Note that for modal logic formulas ϕ we have $\mathcal{M}, \{w\} \models \phi$ iff $\mathcal{M}, w \models \phi$ (where in the latter case, $\models$ is defined as in any textbook for usual modal logic). In fact it is easy to see that without dependence or independence atom, our logic has the so called *flatness property*, stating that team semantics and usual semantics essentially do not make a difference:

Lemma 2.4 *For every* ML*-formula ϕ and all models M and teams T, $M, T \models \phi$ iff $M, w \models \phi$ for all $w \in T$.*

Another simple observation is that the empty team satisfies all formulas.

Lemma 2.5 *For every* MIL*-formula ϕ and all models M, $M, \emptyset \models \phi$.*

Team semantics and independence atoms together will lead to a richer expressive power, as we will prove in Section 6. However, we will also show that over teams T consisting of one world only, ML and MIL have the same expressive power.

Definition 2.6 Formulas φ and φ' are *equivalent on singletons*, if for every model M and every $w \in M$, we have $M, \{w\} \models \varphi$ if and only if $M, \{w\} \models \varphi'$.

Note that on singleton teams, the independence atom trivially always evaluates to true. However, using the modal operators $\Box$ and $\Diamond$, a formula that is evaluated on singleton teams as a starting point clearly is able to talk about nontrivial teams as well. An example for this is the formula constructed in Section 5: The formula is evaluated on a singleton team—the starting point of the protocol—but specifies independence properties for much larger teams, namely subsets of all possible protocol outcomes.

3 Complexity Results

In this section we will study the computational complexity of the model checking and the satisfiability problem for MIL. In [8], it was observed that in first-order team semantics, $=(\vec{p}, q)$ is equivalent to $q \perp_{\vec{p}} q$. This observation clearly carries over to MIL, and hence in particular shows that MIL is a generalization of MDL.

Lemma 3.1 *Let $\mathcal{M}$ be a model and T a team over $\mathcal{M}$, let $\vec{p}$ and q be variables. Then $\mathcal{M}, T \models =(\vec{p}, q)$ if and only if $\mathcal{M}, T \models q \perp_{\vec{p}} q$.*

Proof. Assume that $\mathcal{M}, T \models q \perp_{\vec{p}} q$. We show that $\mathcal{M}, T \models =(\vec{p}, q)$ holds. Let $w, w' \in T$ be such that they agree on the values of $\vec{p}$. We need to show that

$$w \in \pi(q) \Leftrightarrow w' \in \pi(q).$$

Suppose, for a contradiction, that w and w' do not agree on q. By the assumption that $q \perp_{\vec{p}} q$ is satisfied, there exists a world w'' which agrees on $\vec{p}$ with w and w' and is such that $w'' \in \pi(q) \Leftrightarrow w \in \pi(q)$ and $w'' \in \pi(q) \Leftrightarrow w' \in \pi(q)$ which is a contradiction.

The converse implication is proved analogously. □

We now define the two decision problems whose complexity we wish to study, namely the model checking and the satisfiability problem for modal independence logic.

Problem:	MIL-SAT
Input:	MIL formula ϕ
Question:	Does there exists a Kripke model $\mathcal{M}$ and a team $T \neq \emptyset$ with $\mathcal{M}, T \models \phi$?

Problem:	MIL-MC
Input:	Kripke model $\mathcal{M}$, team T and MIL formula ϕ
Question:	$\mathcal{M}, T \models \phi$?

The corresponding problems for modal dependence logic are denoted by MDL-SAT and MDL-MC.

It is easy to see that model checking for MIL is not more difficult than model checking for MDL, namely NP-complete.

Theorem 3.2 MIL-MC *is* NP*-complete.*

Proof. The lower bound follows immediately from Lemma 3.1 and NP-completeness of MDL-MC [6]. The upper bound follows from a simple extension of the well-known model checking algorithm for modal logic, see Algorithm 1.□

Algorithm 1 NP algorithm for MIL-MC

function MILMC$(\mathcal{M}, T, \phi)$
  if $\phi = \Box\psi$ **then**
    return MILMC$(M, R(T), \psi)$
  if $\phi = \Diamond\psi$ **then**
    existentially guess $T' \in R\langle T\rangle$
    return MILMC(M, T', ψ)
  else if $\phi = \psi_1 \wedge \psi_2$ **then**
    return MILMC(M, T, ψ_1) **and** MILMC(M, T, ψ_2)
  else if $\phi = \psi_1 \vee \psi_2$ **then**
    existentially guess $T_1 \cup T_2 = T$
    return MILMC(M, T_1, ψ_1) **and** MILMC(M, T_2, ψ_2)
  else if $\phi = p$ **then**
    for $s \in T$ **do**
      if $p \notin \pi(s)$ **then**
        return false
    return true
  else if $\phi = \overline{p}$ **then**
    for $s \in T$ **do**
      if $p \in \pi(s)$ **then**
        return false
    return true
  else if $\phi = \vec{p} \perp_{\vec{r}} \vec{q}$ **then**
    for $s \in T$ **do**
      for $s' \in T$ **do**
        if $\pi(s) \cap \vec{r} = \pi(s') \cap \vec{r}$ **then**
          found $\leftarrow$ **false**
          for $s'' \in T$ **do**
            agreeP $\leftarrow \pi(s'') \cap \vec{p} = \pi(s) \cap \vec{p}$
            agreeQ $\leftarrow \pi(s'') \cap \vec{q} = \pi(s') \cap \vec{q}$
            agreeR $\leftarrow \pi(s'') \cap \vec{r} = \pi(s) \cap \vec{r}$
            if agreeP **and** agreeQ **and** agreeR **then**
              found $\leftarrow$ **true**
          if not found **then**
            return false
    return true

Next we will consider the complexity of the satisfiability problem MIL-SAT for modal independence logic. From Lemma 3.1 and the hardness of MDL-SAT

for nondeterministic exponential time [15] we immediately obtain the following lower bound:

Lemma 3.3 MIL-SAT *is* NEXP*-hard.*

In order to show containment in NEXP, we need to recall the following classical result. Recall that the so-called Gödel–Kalmár–Schütte prefix class $[\exists^*\forall^2\exists^*, all]$ contains sentences of FO, in a relational vocabulary without equality, which are in prenex normal form and have a quantifier prefix of the form $\exists^*\forall^2\exists^*$.

Proposition 3.4 ([3]) *Satisfiability of formulas in prefix class* $[\exists^*\forall^2\exists^*, all]$ *can be decided in* $\mathrm{NTIME}(2^{O(n/\log n)})$.

Next we will show that MIL-SAT $\in$ NEXP with the help of Proposition 3.4. We will first define a variant of the standard translation of ML into FO that maps MIL-formulas to formulas of monadic existential second-order logic. For a Kripke structure (W, R, π), and a team $T \subseteq W$, we denote by $(W, \{A_i\}_i, R, T)$ the first-order structure of vocabulary $\{R, T\} \cup \{A_i\}_{i\in\mathrm{N}}$ encoding (W, R, π), where relation symbols R and T are interpreted by the accessibility relation R and the team T, and A_i is interpreted by the set $\{w \in W \mid p_i \in \pi(w)\}$.

Lemma 3.5 *For any formula* $\phi \in$ MIL *there is a sentence* ϕ^* *of monadic existential second-order logic of the form*

$$\exists Y_1 \dots \exists Y_m \forall x \forall y \exists z_1 \dots \exists z_k \theta, \tag{1}$$

where θ *is quantifier-free, and such that for all* (W, R, π) *and* T *it holds*

$$(W, R, \pi), T \models \phi \Leftrightarrow (W, \{A_i\}_{i\in\mathrm{N}}, R, T) \models \phi^*.$$

Proof. We first define an auxiliary translation $\phi \mapsto \phi'$ for which correctness is obvious and then indicate how to go from ϕ' to ϕ^*.

(i) Suppose ϕ is p_i. Then ϕ' is defined as

$$\phi' := \forall x(T(x) \to A_i(x)).$$

(ii) Suppose ϕ is $\overline{p}_i$. Then ϕ' is defined as

$$\phi' := \forall x(T(x) \to \neg A_i(x)).$$

(iii) Suppose ϕ is $\psi_1 \vee \psi_2$. Then ϕ' is defined as

$$\phi' := \exists Y_1 \exists Y_2 (\forall x(T(x) \leftrightarrow (Y_1(x) \vee Y_2(x))) \wedge \psi_1'(T/Y_1) \wedge \psi_2'(T/Y_2)).$$

(iv) Suppose ϕ is $\psi_1 \wedge \psi_2$. Then ϕ' is defined as

$$\phi' := \psi_1' \wedge \psi_2'.$$

(v) Suppose ϕ is $\Diamond\psi$. Then ϕ' is defined as

$$\phi' := \exists Y (\forall x(T(x) \to \exists z(Y(z) \wedge E(x, z)) \wedge (Y(x) \to \exists u(T(u) \wedge E(u, x)))) \wedge \psi'(T/Y)).$$

(vi) Suppose ϕ is $\Box\psi$. Then ϕ' is defined as

$$\phi' := \exists Y(\forall x\forall y(((T(x) \wedge E(x,y)) \to Y(y)) \wedge (Y(x) \to \exists z(T(z) \wedge E(z,x)))) \wedge \psi'(T/Y)).$$

(vii) Suppose ϕ is $\vec{p}_1 \perp_{\vec{p}_2} \vec{p}_3$. Then ϕ' is defined as

$$\phi' := \forall x\forall y((T(x) \wedge T(y) \wedge EQ_{\vec{p}_2}(x,y)) \to \exists z(T(z) \wedge EQ_{\vec{p}_2}(x,z) \wedge EQ_{\vec{p}_1}(x,z) \wedge EQ_{\vec{p}_3}(y,z))),$$

where $EQ_{\vec{p}_i}(v,w)$ is a shorthand for the formula

$$\bigwedge_{p_j \in \vec{p}_i} A_j(v) \leftrightarrow A_j(w).$$

It remains to define the translation $\phi \mapsto \phi^*$. This translation is defined by modifying the above clauses by essentially moving all quantifiers to the left of the formula, and by possibly renaming some of the bound variables. We will indicate these modifications by considering the case of disjunction. The other cases are analogous. Assume that ψ_1^* and ψ_2^* are defined already:

$$\psi_i^* = \exists \bar{Y}_i \forall x\forall y\exists \vec{z}_i \theta_i,$$

where θ_i is quantifier free, and $\psi_i^* \equiv \psi_i'$. By renaming of bound variables, we may assume that $\bar{Y}_2 = Y_3 \dots Y_k$, and $\bar{Y}_1 = Y_{k+1} \dots Y_m$, and that $\vec{z}_1$ and $\vec{z}_2$ do not have any common variables either. Then $(\psi_1 \vee \psi_2)^*$ is defined by replacing ψ_i' by ψ_i^* in the definition of $(\psi_1 \vee \psi_2)'$ (see clause iii), and by extending the scopes of the quantifiers:

$$(\psi_1 \vee \psi_2)^* := \exists Y_1 \dots \exists Y_m \forall x\forall y\exists \vec{z}_2\exists \vec{z}_1((T(x) \leftrightarrow (Y_1(x) \vee Y_2(x))) \wedge \theta_1(T/Y_1) \wedge \theta_2(T/Y_2)).$$

□

Theorem 3.6 MIL-SAT *is in* NEXP.

Proof. Let $\phi \in$ MIL. Then ϕ is satisfiable by a Kripke model $\mathcal{M}$ and a team $T \neq \emptyset$ if and only if $\phi^* \wedge \exists w T(w)$ is satisfiable. This follows from the previous lemma and the fact that there is a 1-1 correspondence with Kripke structures (W, R, π), and teams T for ϕ and $\{R, T\} \cup \{A_i\}_{1 \leq i \leq n}$-structures $(W, \{A_i\}_{1 \leq i \leq n}, R, T)$, where n is large enough such that all p_i appearing in ϕ satisfy $i \leq n$.

Recall now that ϕ^* has the form

$$\exists Y_1 \dots \exists Y_m \forall x\forall y\exists z_1 \dots \exists z_k \theta,$$

hence $\phi^* \wedge \exists w T(w)$ is logically equivalent to

$$\exists Y_1 \dots \exists Y_m \forall x\forall y\exists z_1 \dots \exists z_k \exists w(\theta \wedge T(w)),$$

which is satisfiable if and only if the first-order sentence

$$\forall x\forall y\exists z_1 \dots \exists z_k \exists w(\theta \wedge T(w)) \tag{2}$$

of vocabulary $\{Y_1, \ldots, Y_m\} \cup \{R, T\} \cup \{A_i\}_{1 \leq i \leq n}$ is satisfiable. The sentence (2) is contained in prefix class $[\exists^* \forall^2 \exists^*, all]$, hence the satisfiability of it, and also of $\phi^* \wedge \exists w T(w)$, can be decided in time $\mathrm{NTIME}(2^{O(|\phi^*|)})$. The claim now follows from the fact the mapping $\phi \mapsto \phi^*$ can be computed in time polynomial in $|\phi|$. □

Corollary 3.7 MIL-SAT *is* NEXP*-complete.*

It is interesting to note that Theorem 3.6 and Lemma 3.1 directly imply the result of Sevenster [15] that MDL-SAT is contained in NEXP. On the other hand, it seems that the original argument of Sevenster does not immediately generalize to MIL.

Corollary 3.8 MDL-SAT *is* NEXP*-complete.*

4 Generalized Dependency Notions

MIL can be seen as an extension of modal logic with team semantics by the independence atom—let us denote such an extension by $\mathsf{ML}(\perp)$. Similarly, we can extend modal logic with other atoms, so-called *generalized dependence atoms*, which we define now.

Definition 4.1 Let $\mathcal{M} = (W, R, \pi)$ be a Kripke model and $T = (w_1, \ldots, w_m)$ be an arbitrary sequence of elements of W, i.e. essentially an ordered team over $\mathcal{M}$; in the following we will simply refer to these as *teams* when no confusion may arise. Then for any propositional variable p, $T(p)$ is defined as the tuple $(s_1, \ldots, s_m)$, where s_i for $1 \leq i \leq m$ is defined as:

$$s_i = \begin{cases} 1 & w_i \in \pi(p) \\ 0, & \text{otherwise.} \end{cases}$$

For a set of propositions $\vec{q} = (q_1, \ldots, q_k)$, we define $T(\vec{q})$ analogously as $(T(q_1), \ldots, T(q_k))$.

Similar to Kuusisto's [13] definition of generalized first-order dependence atoms we give a definition of generalized modal dependence atoms. In the following, a set of matrices D is invariant under permutations of rows, if for every matrix $M \in D$, if M' is obtained from M by permuting M's rows, then M' is an element of D as well.

Definition 4.2 Let D be a set of Boolean n-column matrices that is invariant under permutation of rows. The semantics of the *generalized dependence atom defined by* D is given as follows:

Let $\mathcal{M}$ be a Kripke model, T be a team over $\mathcal{M}$ and $p_1, \ldots, p_n$ atomic propositions. Then

$$\mathcal{M}, T \models D(p_1, \ldots, p_n) \iff \langle T(p_1), \ldots, T(p_n) \rangle \in D.$$

The *width* of D is defined to be n.

Note that for simplicity we do not distinguish in notation between the logical atom D and the set D of Boolean matrices.

The Boolean matrix $\langle T(p_1), \ldots, T(p_n)\rangle$ contains one column for each of the variables $p_1, \ldots, p_n$; each row of the matrix corresponds to one world from T. The entry for variable p_i and world $w \in T$ is 1 if and only if the variable p_i is satisfied in the world w. We require that D is invariant under permutation of rows in order to ensure that whether $\mathcal{M}, T \models D(p_1, \ldots, p_n)$ holds does not depend on the ordering of the worlds in T that is used in computing the tuple $T(p)$.

In the following we will mainly be interested in generalized dependence atoms *definable* by first-order formulae. For this purpose let D be an atom of width n as above, and ϕ be a first-order sentence over signature $\langle A_1, \ldots, A_n\rangle$. Then ϕ *defines* D if for all Kripke models $\mathcal{M} = (W, R, \pi)$ and teams T over $\mathcal{M}$,

$$\mathcal{M}, T \models D(p_1, \ldots, p_n) \iff \mathcal{A} \models \phi,$$

where $\mathcal{A}$ is the first-order structure with universe T and relations $A_i^{\mathcal{A}}$ for $1 \leq i \leq n$, where for all $w \in T$, $w \in A_i^{\mathcal{A}} \Leftrightarrow p_i \in \pi(w)$.

We say that a generalized dependence atom D is FO-*definable* if there exists a FO-formula ϕ defining D as above. Strictly speaking, the dependence atoms considered in the literature are *families* of dependence atoms for different width, e. g., the simple dependence $=(p_1, \ldots, p_n)$ is defined for arbitrary values of n. Let us say that such a family is (P-uniformly) FO-definable if there exists a family of defining first-order formulae ϕ_n such that ϕ_n defines the atom of width n and the mapping $1^n \mapsto \langle \phi_n \rangle$ is computable in polynomial time; that is, an encoding of formula ϕ_n is computable in time polynomial in n. Note that in particular this implies that $|\phi_n| = p(n)$ for some polynomial p.

As examples let us show how to define some well-studied generalized dependence atoms as follows.

$$\begin{array}{ll} =(\vec{p}, q) & \Leftrightarrow \forall w \forall w' \ ((\bigwedge_{1\leq i\leq n} A_{p_i}(w) \leftrightarrow A_{p_i}(w')) \rightarrow (A_q(w) \leftrightarrow A_q(w'))) \\ \vec{p} \subseteq \vec{q} & \Leftrightarrow \forall w \exists w' \ (\bigwedge_{1\leq i\leq n} A_{p_i}(w) \leftrightarrow A_{q_i}(w')) \\ \vec{p} \mid \vec{q} & \Leftrightarrow \forall w \forall w' \ (\bigvee_{1\leq i\leq n} A_{p_i}(w) \leftrightarrow \neg A_{q_i}(w')) \end{array}$$

The latter two so-called *inclusion* and *exclusion* atoms were introduced by Galliani in [7]. In particular all above atoms are FO-definable. The independence atom $\vec{p}_1 \perp_{\vec{q}} \vec{p}_2$ is also FO-definable in the obvious way. [1]

We use $\mathsf{ML}(D)$ to denote the extension of ML by a generalized dependence atom D. We will next show that our complexity upper bounds from Section 3 can be generalized to cover $\mathsf{ML}(D)$ for certain FO-definable dependence atoms D. For model checking, we simply use the fact that first-oder formulas can be verified in polynomial time and obtain the following corollary:

[1] Since the FO-formula ϕ may only depend on the width, we restrict ourselves to occurrences of $\vec{p}_1 \perp_{\vec{q}} \vec{p}_2$ where $|\vec{p}_1| = |\vec{p}_2| = |\vec{q}|$, if these sets are nonempty, which we can always assume without loss of generality by repeating variable occurrences, the case that one of these sets is empty can then be encoded into widths that are not multiples of 3 in a straightforward manner.

Corollary 4.3 *Let D be a P-uniformly* FO*-definable generalized dependence atom. Then* $\mathsf{ML}(D)$-MC *is in* NP*.*

For the satisfiability problem, we generalize the proof of Theorem 3.5 in case (vii) to dependence atoms which are definable by a $[\exists^*\forall^2\exists^*]$ formula.

Corollary 4.4 *Let D be a generalized dependence atom that is P-uniformly* FO*-definable by a (family of) first-order formula(e) in the prefix class* $[\exists^*\forall^2\exists^*]$*. Then* $\mathsf{ML}(D)$-SAT *is in* NEXP*.*

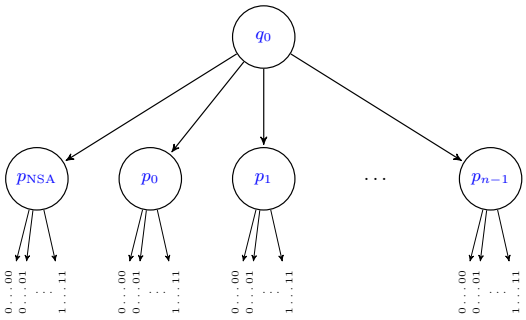
We remark that the prefix class $[\exists^*\forall^2\exists^*]$ is expressive enough to define to the best of our knowledge all generalized dependence atoms studied in the literature so far.

5 Example: The Dining Cryptographers

The dining cryptographers [4], a standard example for anonymous broadcast, is the following problem: A group of cryptographers $\{c_0, \ldots, c_{n-1}\}$ with $n \geq 3$ sit in a restaurant, where c_i sits between c_{i-1} and c_{i+1}. (Indices of the cryptographers are always modulo n, and i always ranges over $0, \ldots, n-1$). After dinner, it turns out that someone already paid. There are only two possibilities: Either one of the cryptographers secretly paid, or the NSA did. Naturally, they want to know which of these is the case, but without revealing the paying cryptographer if one of them paid. They use the following protocol (in the protocol, $\oplus$ defines the exclusive-or of two bits, where $b_1 \oplus b_2 = b_1 + b_2 \bmod 2$):

- For each i, let p_i be 1 iff c_i paid. Each c_i knows the value of p_i, but not of p_j for $j \neq i$. There is at most one i with $p_i = 1$. The protocol computes the value $p_0 \oplus p_1 \oplus \cdots \oplus p_{n-1}$, which is the same as $p_0 \vee p_1 \vee \cdots \vee p_{n-1}$.
- Each adjacent pair $\{c_i, c_{i+1}\}$ computes a random bit $\mathit{bit}_{\{i,i+1\}}$.
- Each c_i publicly announces the value $\mathit{announce}_i = p_i \oplus \mathit{bit}_{\{i,i-1\}} \oplus \mathit{bit}_{\{i,i+1\}}$.
- Then, $p_0 \oplus p_1 \oplus \cdots \oplus p_{n-1} = \mathit{announce}_0 \oplus \mathit{announce}_1 \oplus \cdots \oplus \mathit{announce}_{n-1}$.

The protocol clearly computes the correct answer, the interesting aspect is the *anonymity requirement*: No cryptographer c_i should learn anything about the values p_j for $j \neq i$ except for what follows from the values p_i or the result (if c_i or the NSA paid then c_j did not). The protocol models anonymous broadcast, since the message "1" is, if sent, received by all cryptographers, but the sender remains anonymous. We formalize this using modal independence logic. We start by capturing the protocol in the following Kripke model:



The protocol starts in q_0, the model then branches into states p_{NSA}, p_0, …, p_{n-1}, depending on whether the NSA or some c_i paid. Each of these states has 2^n successor states, for the 2^n possible random bit values, these states are *final*. The relation R is as indicated. We use the following variables:

- p_{NSA} and p_i are true if the NSA, resp. cryptographer c_i paid, i.e., in the states denoted with the same name as the variable and in their successors.
- each of the n variables $bit_{\{i,i+1\}}$ is true in the states where the bit shared between c_i and c_{i+1} is 1.
- each $announce_i$ is true in all final states which satisfy $p_i \oplus bit_{\{i,i-1\}} \oplus bit_{\{i,i+1\}}$ (this encodes that the cryptographers follow the protocol).

For each c_i, we define the set $\vec{k}_i$ of the variables whose values c_i knows after the protocol run as $\vec{k}_i := \{p_i, bit_{\{i,i-1\}}, bit_{\{i,i+1\}}\} \cup \{announce_j \mid j \neq i\}$. Clearly, c_i also knows the value $announce_i$, but since this can be computed from p_i, $bit_{\{i,i-1\}}$ and $bit_{\{i,i+1\}}$, we omit it from $\vec{k}_i$.

The formula expressing the anonymity requirement consists of several parts, one *global* part and then, for each combination of cryptographers, a *local* part. We start with the global part, which merely expresses that none of the individual bits that some cryptographer knows determines the value of any p_i on its own, with the exception that if $p_i = 1$, then of course cryptographer c_i knows that $p_j = 0$ for all $j \neq i$. The global part φ_g is as follows:

$$\varphi_g = \bigwedge_{\substack{i\in\{0,\dots,n-1\},\\ v\in\{bit_{\{i,i+1\}},announce_i\},\\ k\in\{0,\dots,n-1\}}} \Diamond\Diamond(v \wedge p_k) \wedge \Diamond\Diamond(v \wedge \overline{p_k}) \wedge \Diamond\Diamond(\overline{v} \wedge p_k) \wedge \Diamond\Diamond(\overline{v} \wedge \overline{p_k}) \\ \wedge \textstyle\bigwedge_{i\neq j} \Diamond\Diamond(p_i \wedge \overline{p_j}) \wedge \Diamond\Diamond(\overline{p_i} \wedge p_j) \wedge \Diamond\Diamond(\overline{p_i} \wedge \overline{p_j})$$

The first line of the formula requires that, for every variable v of the $bit_{\{i,i+1\}}$ or $announce_i$-variables, and every cryptographer c_k, every combination of truth values of v and p_k appears. This encodes that the value of a *single* variable v does not give away any information about the value of any p_k. The second line is a similar requirement for the value p_i: If c_i paid, then she knows that c_j did not pay, for $i \neq j$. However, the combination "$p_i \wedge p_j$" for $i \neq j$ should be the only one not appearing. Hence the formula requires that

all other combinations appear in some final state. The global part φ_g hence ensures that each individual bit that c_i knows does not tell him whether c_j paid, unless of course $i = j$ or $p_i = 1$.

The more interesting part is to encode that even the *combination* of the above bits does not lead to additional knowledge; this is where the independence atom is crucial. We introduce some notation to enumerate the variables in $\vec{k}_i$:

- for each i, let $\vec{k}_i = \{v_1^i, \ldots, v_{n+2}^i\}$ be an arbitrary enumeration of the variables in $\vec{k}_i$,
- for $j \leq k$, let $V_{j\to k}^i = \{v_j^i, \ldots, v_k^i\}$,
- let $V_j^i = V_{j\to j}^i$.

We now use modal independence logic to express that if each single variable from $\vec{k}_i$ does not tell c_i anything about the value of p_k, then their combination does not, either. This is achieved with the following formula:

$$\varphi^{i,k} = \Box\Box\left((V_1^i \perp_{p_k} V_2^i) \wedge (V_{1\to 2}^i \perp_{p_k} V_3^i) \wedge \cdots \wedge (V_{1\to n+1}^i \perp_{p_k} V_{n+2}^i)\right).$$

This formula requires that for each j, each pair of variable assignments I_1 to $V_{1\to j-1}^i$ and I_2 to V_j^i that is "locally compatible" with some truth value $P(p_k)$—in other words, neither of these assignments by itself implies that the actual value of p_k is not $P(p_k)$—is also compatible with that value for the combination of I_1 and I_2, i.e., there is some state satisfying $I_1 \cup I_2 \cup P$ (where the notion of a state satisfying a propositional assignment is defined as expected and the union of these assignments is well-defined since their domains are disjoint). As a consequence, the formula requires that for each $I\colon \vec{k}_i \to \{0,1\}$ and each $P\colon \{p_k\} \to \{0,1\}$, if for each $v \in \vec{k}_i$, there is a world w such that $w \models I\big|_{\{v\}}$ and $w \models P$, then there is a world w such that $w \models I \cup P$.

The following proposition formally states that our above-developed formulas indeed express the anonymity property of the protocol as intended. From a single cryptographer c_i's point of view, it says that every observation I which can arise when c_i follows the protocol, as long as some cryptographer different from c_i paid for the dinner, then for every k different from i, both possibilities—c_k paid for the dinner, or c_k did not pay—cannot be ruled out by the observation I. We say that an assignment $I\colon \vec{k}_i \cup \{announce_i\} \to \{0,1\}$ is *consistent* if i follows the protocol, i.e., if $I(announce_i) = I(p_i) \oplus I(bit_{\{i,i-1\}}) \oplus I(bit_{\{i,i+1\}})$. Note that in the models we are interested in, only consistent assignments appear in final states.

Proposition 5.1 *If a Kripke model $M = (W, R, \pi)$ satisfies the formula $\varphi_g \wedge \bigwedge_{i,k\in\{0,\ldots,n-1\}, i\neq k} \varphi^{i,k}$ at the world q_0, then the team $T = R(R(\{q_0\}))$ satisfies the following condition: For each $i \neq k \in \{0,\ldots,n-1\}$ and each consistent $I\colon \vec{k}_i \cup \{announce_i\} \to \{0,1\}$ with $I(p_i) = 0$ and $\oplus_{j=0}^{n-1} I(announce_j) = 1$, there are worlds $w_1^I, w_2^I \in T$ with $w_1^I \models I, p_k$ and $w_2^I \models I, \overline{p_k}$.*

We omit the easy proof; the proposition immediately follows from the se-

mantics of the independence atom.

Our discussion only treats the anonymity property of the protocol. For a complete treatment, one also has to address other aspects as e.g., correctness, we omit this discussion here.

Note that is is also possible to express the anonymity requirement using epistemic logic (see, e.g., [1,14]). Usually, this is done by explicitly introducing, for each participant a of the protocol, a modality R_a that models her information in the sense that two worlds are related with R_a are indistinguishable for the participant a. (Such a modality is then necessarily an equivalence relation). The main difference to our modelling is that we do not use the modality of the Kripke model to represent knowledge, but to express branching time. In particular, our approach only uses a single modality, and reasons about knowledge of the protocol's principals using the involved propositional variables.

6 Expressiveness

We now compare the expressiveness of MIL and classical modal logic, which we abbreviate with ML. We show that MIL is strictly more expressive than ML on teams (simply because MIL is not downwards closed, i.e., from $\mathcal{M}, T \models \varphi$ and $T' \subseteq T$, it does not follow that $\mathcal{M}, T' \models \varphi$), but that their expressiveness coincides on singleton teams. However, on singletons, MIL is exponentially more succinct than ML. We then study the expressiveness of MIL with a generalized dependence atom as introduced in Section 4 instead of the independence atom.

6.1 Expressiveness of MIL and ML

Clearly, since MIL is not downward-closed, we obtain the following:

Proposition 6.1 *There is an* MIL*-formula* φ_{MIL} *such that there is no* ML*-formula* φ_{ML} *such that the following property holds:* $M, T \models \varphi_{\mathsf{MIL}}$ *if and only if* $M, T \models \varphi_{\mathsf{ML}}$ *for all models* M *and all teams* T*.*

Proof. This is true for every formula φ_{MIL} that is not downwards closed: In this case we have teams $T' \subsetneq T$ of the same model M with $M, T \models \varphi_{\mathsf{MIL}}$ and $M, T' \not\models \varphi_{\mathsf{MIL}}$. However, for any modal formula φ_{ML}, clearly if $M, w \models \varphi_{\mathsf{ML}}$ for all $w \in T$, then the same is true for all $w \in T'$ as $T' \subseteq T$. An easy example for a formula that is not downwards closed is $x \perp_\emptyset y$. This formula is satisfied on a team T in which every combination of truth values of x and y is realized in some world, but not on its subset T' containing only worlds w and w' with assignments $x \wedge y$ and $\overline{x} \wedge \overline{y}$, respectively. □

The proposition remains true when you replace ML by classical modal logic extended with a global modality, or by MDL, since these logics remain downward-closed. In [15], it was shown that MDL is as expressive as classical modal logic on singletons. Therefore, a natural question to ask is whether on singletons, MIL is still more expressive than ML. We show that this is not the case, but we will also see that MIL is exponentially more succinct than ML, even on singletons. For our proof, we use bisimulations, which are a well-established tool to compare expressiveness of different concepts. We recall the classical

definition of bisimulation for modal logic:

Definition 6.2 Let $M = (W, R, \pi)$ and $M' = (W', R', \pi')$ be Kripke models. A relation $Z \subseteq W \times W'$ is a *modal bisimulation* if for every $(w, w') \in Z$, the following holds:

- $\pi(w) = \pi'(w')$, i.e., w and w' satisfy the same propositional variables,
- if u is an R-successor of w, then there is an R'-successor u' of w' such that $(u, u') \in Z$ (forward condition),
- if u' is an R'-successor of w', then there is an R-successor u of w such that $(u, u') \in Z$ (backward condition).

It is well-known and easy to see that modal logic is invariant under bisimulation, i.e., if Z is a bisimulation and $(w, w') \in Z$, then w and w' satisfy the same modal formulas. We now "lift" this property to modal independence logic by considering a bisimulation Z as above on the team level:

Definition 6.3 Let $M = (W, R, \pi)$ and $M' = (W', R', \pi')$ be models, let $T \subseteq W$ and $T' \subseteq W'$ be teams. Let $Z \subseteq W \times W'$ be a modal bisimulation. Then T and T' are Z-bisimilar if the following is true:

- for each $w \in T$, there is a $w' \in T'$ such that $(w, w') \in Z$,
- for each $w' \in T'$, there is a $w \in T$ such that $(w, w') \in Z$.

We now show that on the team level, bisimulation for modal independence logic plays the same role as it does on the world level for modal logic: Simply stated, bisimilar teams satisfy the same formulas. Due to Lemma 3.1, the result also applies to modal dependence logic. This lemma may be of independent interest (for example, it implies a "family-of-trees"-like model property), we use it to compare the expressiveness of MIL and ML.

Lemma 6.4 *Let $M = (W, R, \pi)$ and $M' = (W', R', \pi')$ be Kripke models, let $T \subseteq W$ and $T' \subseteq W'$ be teams that are Z-bisimilar for a modal bisimulation Z. Then for any MIL-formula φ, we have that $M, T \models \varphi$ if and only if $M', T' \models \varphi$.*

Proof. We show the lemma by induction on φ. Clearly it suffices to show that if $M, T \models_{\mathsf{MIL}} \varphi$, then $M', T' \models_{\mathsf{MIL}} \varphi$. Hence assume $M, T \models_{\mathsf{MIL}} \varphi$.

- Let $\varphi = x$ for some propositional variable x, and let $w' \in T'$. Since T and T' are Z-bisimilar, there is a world $w \in T$ with $(w, w') \in Z$. Since $M, T \models_{\mathsf{MIL}} x$, the variable x is true at w in M. Since Z is a modal bisimulation, it follows that x is true at w' in M', and hence every world $w' \in T$ satisfies x. Therefore, it follows that $M', T' \models \varphi$.
- If $\varphi = \neg x$, the proof is the same as above.
- Let $\varphi = \varphi_1 \wedge \varphi_2$. This case trivially follows inductively.
- Let $\varphi = \varphi_1 \vee \varphi_2$. Since $M, T \models \varphi$, it follows that $T = T_1 \cup T_2$ for teams T_1 and T_2 with $M, T_1 \models \varphi_1$ and $M, T_2 \models \varphi_2$. We define teams T_1' and T_2' as follows:
  · $T_1' = \{w' \in T' \mid (w, w') \in Z \text{ for some } w \in T_1\}$,

· $T_2' = \{w' \in T' \mid (w, w') \in Z \text{ for some } w \in T_2\}$.

We prove the following:

(i) $T' = T_1' \cup T_2'$

(ii) T_1 and T_1' are Z-bisimilar,

(iii) T_2 and T_2' are Z-bisimilar.

By induction, it then follows that $M', T_1' \models \varphi_1$ and $M', T_2' \models \varphi_2$, which, since $T' = T_1' \cup T_2'$ implies that $M', T' \models \varphi$. We prove these points:

(i) By construction, $T_1' \cup T_2' \subseteq T'$. Hence let $w' \in T'$. Since T and T' are Z-bisimilar, there is some $w \in T$ such that $(w, w') \in Z$. Since $T = T_1 \cup T_2$, we can, without loss of generality, assume that $w \in T_1$. By definition of T_1', it follows that $w' \in T_1'$.

(ii) First let $w \in T_1 \subseteq T$. Since T and T' are Z-bisimilar, there is some $w' \in T'$ such that $(w, w') \in Z$. Due to the definition of T_1', it follows that $w' \in T_1'$. For the converse, assume that $w' \in T_1'$. By definition, there is some $w \in T_1$ such that $(w, w') \in Z$.

(iii) This follows with the same proof as for T_1 and T_1'.

Hence $M', T' \models \varphi$ as required.

- Let $\varphi = \Diamond\psi$. Since $M, T \models \varphi$, there exists a team $U \subseteq R(T)$ such that for each $w \in T$, the set $u(w) := R(\{w\}) \cap U$ is not empty, and $M, U \models \psi$. We define a corresponding team U' of M' as follows: Start with $U' = \emptyset$ and then for each $(w, w') \in (T \times T') \cap Z$, do the following:

  · For each R-successor v of w that is an element of U, since Z is a modal bisimulation and $(w, w') \in Z$, there is at least one R'-successor v' of w' with $(v, v') \in Z$. Add all such v' to the set U'.

  By construction, U' only contains worlds that are R'-successors of worlds in T'. Hence to show that $M', T' \models \Diamond\psi$, it remains to show that

  (i) for each $w' \in T'$, the team U' contains a world v' that is an R'-successor of w',

  (ii) the teams U and U' are Z-bisimilar.

  The claim then follows by induction, since $M, U \models \psi$. We now show the above two points:

  (i) Let $w' \in T'$. Since T and T' are Z-bisimilar, there is some $w \in T$ with $(w, w') \in Z$. Due to the choice of U, there is some $v \in U$ which is an R-successor of w. By construction of the set U', a world v' that is an R'-successor of w' has been added to U'.

  (ii) By construction, for every R-successor v of some w in T, at least one v' has been added to U' with $(v, v') \in Z$. For the converse, by construction of U', for every v' added to U' there is a $v \in U$ with $(v, v') \in Z$.

  Hence $M', T' \models \Diamond\psi$ as required.

- Now assume that $\varphi = \Box\psi$, and let U be the set of all R-successors of worlds in T, let U' be the set of all R'-successors of worlds in T'. By induction, it suffices to show that U and U' are Z-bisimilar. Hence let $v \in U$ be the R-successor of some $w \in T$. Since T and T' are Z-bisimilar, there is some $w' \in T'$ such that $(w, w') \in Z$. Since v is an R-successor of w, and since Z

is a modal bisimulation, there is some v' which is an R'-successor of w' such that $(v, v') \in Z$. Since v' is an R'-successor of w', it follows that $v' \in U'$. The converse direction follows analogously.

- Let $\varphi = \vec{p_1} \perp_{\vec{q}} \vec{p_2}$. To show that $M', T' \models \varphi$, let $u, u' \in T'$ with the same truth values of the variables in q. Since T and T' are bisimilar, there are worlds $w, w' \in T$ such that $(w, u) \in Z$ and $(w', u') \in Z$. Since Z is a modal bisimulation, the Z-related worlds have the same propositional truth assignment. In particular, w and w' agree on the values for the variables in q. Since $M, T \models \varphi$, there is some world $w'' \in T$ such that
  · $w'' \equiv_{\vec{q}} w' \equiv_{\vec{q}} w$, and since Z is a modal bisimulation it follows that w'' and both u and u' have the same $\vec{q}$-assignment,
  · $w'' \equiv_{\vec{p_1}} w$, and hence w'' and u have the same $\vec{p_1}$-assignment,
  · $w'' \equiv_{\vec{p_2}} w'$, and hence w'' and u' have the same $\vec{p_2}$-assignment.
  Since T and T' are Z-bisimilar, there is a world $u'' \in T'$ such that $(w'', u'') \in Z$. Since Z is a modal bisimulation, w'' and u'' satisfy the same propositional variables, and hence for u'' we have that
  · $u'' \equiv_{\vec{q}} u$, $u'' \equiv_{\vec{q}} u'$
  · $u'' \equiv_{\vec{p_1}} u$
  · $u'' \equiv_{\vec{p_2}} u'$.
  Therefore, $M', T' \models \varphi$ as required.

□

With Lemma 6.4 and an application of van Benthem's Theorem [2], it follows directly that MIL and ML are in fact equivalent in expressiveness *on singletons*. In particular, this implies that the formula we constructed in Section 5 can be rewritten as an equivalent formula of classical modal logic-recall that we only evaluated the formula in a singleton world. We will see shortly, however, that our modelling gives us a *succinctness* advantage over using classical modal logic.

The proof of the following theorem relies on van Benthem's Theorem, which states that ML is expressively equivalent with the bisimulation-invariant fragment of first-order logic.

Theorem 6.5 *For each* MIL*-formula* φ_{MIL}*, there is an* ML*-formula* φ_{ML} *such that* φ_{ML} *and* φ_{MIL} *are equivalent on singletons.*

Proof. Due to Lemma 6.4, we know that MIL is invariant under bisimulation of teams. Since for singleton teams, bisimulation on teams and bisimulation on worlds coincide, it follows that MIL on singletons is invariant under modal bisimulation. Clearly, when evaluating an MIL-formula φ_{MIL} on a singleton team $\{w\}$, all worlds in the model that have a distance from w which exceeds the modal depth (i.e., maximal nesting degree of modal operators) of φ, are irrelevant for the question whether $M, \{w\} \models \varphi$. Therefore, φ_{MIL}, evaluated on singletons, captures a property of Kripke models that is invariant under modal bisimulation and only depends on the worlds that can be reached in at most $md(\varphi_{\mathsf{MIL}})$ steps. Due to [2], such a property can be encoded by a standard modal logic formula φ_{ML}. (The formula φ_{ML} can be obtained, for

example, as the disjunction of formulas $\varphi_{M,w}$ which, for each model M and world w with $M, \{w\} \models_{\mathsf{MIL}} \varphi_{\mathsf{MIL}}$, encodes the finite tree unfolding of M, w up to depth $md(\varphi_{\mathsf{MIL}})$, up to bisimulation, and only taking into account the variables appearing in φ_{MIL}. This unfolding is finite and hence first-order definable, therefore we can apply the result from [2].) □

Since the application of van Benthem's Theorem yields a potentially very large formula, the above result does not give an "efficient" translation from MIL to ML. It turns out that one cannot do much better: MIL is exponentially more succinct than ML.

Theorem 6.6 *There is a family of* MIL*-formulas* $(\varphi_i)_{i\in\mathbb{N}}$ *such that the length of* φ_i *is quadratic in* i*, and for any family of* ML*-formulas* $(\psi_i)_{i\in\mathbb{N}}$ *such that for all* i*,* φ_i *and* ψ_i *are equivalent on singletons, the length of* ψ_i *grows exponentially in* i*.*

Proof. Let φ_i be the formula describing the security property of the dining cryptographers protocol for i cryptographers, as constructed in Section 5, with every sequence $\Diamond\Diamond$ replaced with $\Diamond$. As argued in that section, if φ_i is satisfied at $M, \{w\}$, then the number of propositional assignments appearing in the set of worlds that can be reached from w in one step (two steps for the original formula) is exponential in i. Now let $(\psi_i)_{i\in\mathbb{N}}$ be a family of ML-formulas such that φ_i and ψ_i are locally equivalent for each i. Since ψ_i is locally equivalent to φ_i, we can without loss of generality assume that $md(\psi_i) = 1$ for all i (if ψ_i contains deeper nestings of modal operators, the formula can be simplified since the truth value of ψ_i cannot depend on worlds reachable in 2 or more steps). Therefore, modal operators do not appear nested in ψ_i. Without loss of generality, we can assume that only $\Diamond$ appears in ψ. It is clear that if $M, w \models \psi_i$, then there is a submodel of M which contains w, and in which the number of successors of w is bounded by the number of $\Diamond$-operators appearing in ψ_i. Therefore, ψ_i must have an exponential number of $\Diamond$-operators, which proves the theorem. □

As far as we know, the analogue of Theorem 6.6 for MDL has not yet been showed. Sevenster [15] has shown the analogue of Theorem 6.5 for MDL, but he did not show that going from MDL to ML (over singletons) inevitably leads to an exponential blow-up in the formula size. On the other hand, a closely related result showing that any formula of $\mathsf{ML}(\veebar)$, where $\veebar$ is the classical disjunction) that is logically equivalent to

$$=(p_1, \ldots, p_n, q),$$

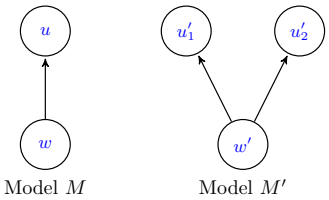
has to have length exponential in n was obtained in [10].

6.2 Expressiveness of ML with Generalized Dependence Atoms

Theorem 6.5 applies not only to MIL, but (with the same proof) to all extensions of MIL with a dependence operator that, evaluated on a team T, depends only on the set of propositional assignments that occur in some world of the team.

This is because the set of assignments is clearly invariant under bisimulation. As examples, operators like those from exclusion logic or inclusion logic can be added to MIL without increasing its expressiveness. Hence, in light of Section 4, a natural question to ask is the following: For which generalized dependence atoms D is the logic $\mathsf{ML}(D)$ as expressive as ML on singletons, and which atoms do in fact add expressiveness?

It is easy to see that there are generalized dependence atoms D that, even on singletons, add expressiveness beyond classical modal logic (and hence, beyond MIL). This is because with no restriction on the dependence operator D, one can express properties that depend on the *number* of worlds in a team, which clearly cannot be done in ML. As an example, consider the following:



Example 6.7 Let D be the relation $\{(0)\}$, and consider the models M and M' above, where the single variable x is false in every world of both models. It is easy to see that $M, \{w\} \models \Box D(x)$, while $M', \{w'\} \not\models \Box D(x)$: For the model M, the set of successors of w is the team $T = \{u\}$, since x is false in u, it follows that $T(x) = (0)$, hence $T \models D(x)$. On the other hand, the set of successors of w' in M' is $T' = \{u_1, u_2\}$, hence $T'(x) = (0, 0)$ and $T' \not\models D(x)$.

Clearly, no ML-formula can distinguish M, w and M', w', since the relation $Z = \{(w, w'), (u, u_1), (u, u_2)\}$ is a bisimulation.

However, as mentioned earlier, the proof of Theorem 6.5 can be generalized to handle the generalized dependence atoms discussed earlier. In general, we obtain the following result: For a D which is definable using first-order formulas without equality, the expressiveness of $\mathsf{ML}(D)$ coincides with that on ML for singletons. For the proof, we show that $\mathsf{ML}(D)$ remains invariant under bisimulation, and then apply the proof of Theorem 6.5 again. On the other hand, clearly if D is not FO-definable, then D cannot be expressed in modal logic, as modal logic can be translated to first-order logic with the standard translation, and if D cannot be expressed without equality, then D cannot be invariant under bisimulation. Hence we obtain the following theorem:

Theorem 6.8 *Let D be a generalized dependence atom. Then the following statements are equivalent:*

(i) *D can be expressed in first-order logic without equality,*

(ii) *$\mathsf{ML}(D)$ and ML are equally expressive over singletons, i.e., for each $\mathsf{ML}(D)$-formula φ_{MIL}, there is an ML-formula φ_{ML} such that φ_{MIL} and φ_{ML} are equivalent on singletons.*

Proof. We first assume that D is FO-definable. As mentioned above, it suffices

to adapt the proof of Lemma 6.4 to $\mathsf{ML}(D)$. Clearly, in the induction, we only need to cover the case that $\varphi = D(p_1, \ldots, p_n)$ for propositional variables $p_1, \ldots, p_n$. Hence assume that $M, T \models D(p_1, \ldots, p_n)$, we show that $M', T' \models D(p_1, \ldots, p_n)$, where T and T' are Z-bisimilar for a modal bisimulation Z.

Let ϕ be the first-order formula defining D. We prove the claim inductively over the formula, where we only cover the key cases explicitly. Let the free variables of ϕ be $\omega_1, \ldots, \omega_t$. We show that if $w_1, \ldots, w_t \in T$ and $w'_1, \ldots, w'_n \in T'$ such that $(w_i, w'_i) \in Z$ for all i, then $\phi(w_1, \ldots, w_n)$ evaluates to true if and only if $\phi(w'_1, \ldots, w'_n)$ does. If ϕ is quantifier-free, the claim is clear: Since Z is a modal bisimulation, w_i and w'_i satisfy the same propositional variables, and due to the choice of ϕ, the truth value of ϕ only depends on the propositional assignments of the worlds instantiating the variables $\omega_1, \ldots, \omega_n$. The second relevant case is when $\phi = \exists\omega\psi(\omega, \omega_1, \ldots, \omega_t)$. If $M, T \models \phi$, then there is a world $w \in T$ such that $\psi(w, w_1, \ldots, w_t)$ is true. Since T and T' are Z-bisimilar, it follows that there is a world $w' \in T'$ such that $(w, w') \in Z$. Due to induction, it follows that if $\psi(w, w_1, \ldots, w_t)$ is true, then so is $\psi(w', w'_1, \ldots, w'_t)$. This completes the proof that $\mathsf{ML}(D)$ and ML are equally expressive over singletons.

For the converse, assume that $\mathsf{ML}(D)$ is as expressive as ML over singletons. In particular, then for every sequence $x_1, \ldots, x_n$ of variables, there is a modal formula φ such that for every model M and every world $w \in M$, we have that $M, w \models \varphi$ if and only if $M, \{w\} \models \Box D(x_1, \ldots, x_n)$. By the standard translation from modal logic to first-order logic, this implies that $D(x_1, \ldots, x_n)$ can be expressed as a FO-formula ϕ_D.

Since whether $\mathcal{M}, T \models D(x_1, \ldots, x_n)$ is invariant under bisimulation (this can be seen by adding a single world w_0, connected to all $w' \in T$ and evaluating the formula $\Box D(x_1, \ldots, x_n)$ at the singleton team $\{w_0\}$—since $\mathsf{ML}(D)$ on singletons is equivalent to ML, which is invariant under bisimulation, it follows that $\mathsf{ML}(D)$ on singletons is invariant under bisimulation as well), it follows that whether $\mathcal{M}, T \models D(x_1, \ldots, x_n)$ is in particular invariant under adding/removing identical copies of worlds in T. Therefore, ϕ_D can be rewritten into a formula without equality. □

7 Conclusion and Open questions

In this paper we introduced modal independence logic MIL and settled the computational complexity of its satisfiability and model checking problem. Furthermore we compared the expressivity of MIL with that of classical modal logic. It turned out that most of our results can be generalized to modal logic extended with various generalized dependence atoms.

The generalization of bisimilarity defined in this article (Definition 6.3) has been recently studied in [11], where it was shown that the so-called Modal Team Logic (MTL) extending MDL by the classical negation can define exactly the FO-definable bisimulation invariant properties of Kripke structures and teams. It was also observed that MIL is a proper sublogic of MTL. It is an open question whether MIL corresponds to a natural subclass of the FO-definable bisimulation invariant properties.

In this article we pinned down the complexity of the satisfiability and the model checking problems of MIL, but the complexity of the validity problem of MIL is still open. It worth noting that the complexity of this problem for the propositional fragment of MIL has been studied in [9].

Acknowledgement

The first author was supported by the Academy of Finland grants 292767 and 275241.

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